

Application of Information Theory, Lecture 5

Channel Capacity, Isoperimetric inequality

Handout Mode

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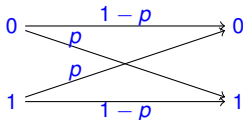
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Part I

Channel Capacity

The problem

- ▶ We want to send a message $\mathbf{x} = (x_1, \dots, x_m) \in \{0, 1\}^m$, but the communication channel is **faulty**
- ▶ Each bit is (independently) flipped w.p. p (e.g., 0.1)



- ▶ (expected) **Error rate** is p
- ▶ Such “channel” is called **Binary Symmetric Channel (BSC)**
- ▶ When sending m bits, we have $\approx pm$ errors
- ▶ Can we send bits with smaller error?

Solution

- ▶ Obvious solution is “error correction codes (ECC)”
- ▶ Most simple example: send each bit three times, and take majority
- ▶ Error happens if the channel errs at least twice
- ▶ For $p = .1$: happens w.p. $3p^2(1 - p) + p^3 = 3 \cdot .01 \cdot .9 + .001 = .028$
- ▶ Error rate: .028
- ▶ **Transmission rate**: $1/3$ (i.e., # of bits recovered / #of bits transmitted)
- ▶ We reduced the error rate by reducing the transmission rate.
- ▶ Can we reduce the error rate, **without** reducing the transmitting rate too much?
- ▶ Before Shannon it was believed that very small error rate requires very small transmission rate.

Shannon's result

- ▶ Shannon showed that you can reduce the error rate towards 0, without reducing the transmission rate towards 0
- ▶ For any $c < C_p$, exists a code with transmission rate c that is correct w.p.
- ▶ Example: for $p = .1$, $C_p > \frac{1}{2}$.

Hence, for sending $\mathbf{x} = (x_1, \dots, x_m)$, one can send $2m$ bits, such that \mathbf{x} is recovered w.p. close to 1

- ▶ More generally, $\forall p \in [0, 1] \exists C_p$ such that for sending $\mathbf{x} \in \{0, 1\}^m$, one can send $\approx m \cdot C_p$ bits, and \mathbf{x} is recovered w.p. close to 1
- ▶ C_p might be 0 (i.e., for $p = \frac{1}{2}$)
- ▶ A revolution in EE and the whole world

Error correction code

- ▶ Message to send $\mathbf{x} = (x_1, \dots, x_m) \in \{0, 1\}^m$
- ▶ Encoding scheme: $f: \{0, 1\}^m \mapsto \{0, 1\}^n$ ($n > m$)
- ▶ Decoding scheme: $g: \{0, 1\}^n \mapsto \{0, 1\}^m$
- ▶ $\frac{m}{n}$ — transmission rate
- ▶ Sender sends $f(\mathbf{x})$ rather than \mathbf{x}
- ▶ Receiver decodes the message by applying g

$$\underbrace{\mathbf{x}}_{m \text{ bits}} \xrightarrow{\text{encoding}} \underbrace{f(\mathbf{x})}_{n \text{ bits}} \xrightarrow{\text{channel}} \underbrace{f(\mathbf{x}) \oplus Z}_{\text{bitwise XOR}} \xrightarrow{\text{decoding}} g(f(\mathbf{x}) \oplus Z)$$

$Z = (Z_1, \dots, Z_n)$ where Z_1, \dots, Z_n iid $\sim (1 - p, p)$ (i.e., over $\{0, 1\}$ with $\Pr[Z_i = 1] = p$)

- ▶ We hope $g(f(\mathbf{x}) \oplus Z) = \mathbf{x}$
- ▶ ECCs are everywhere
- ▶ ECC Vs compression

Shannon's theorem

Theorem 1

$\forall p \quad \exists C_p, \text{ s.t. } \forall \varepsilon > 0 \quad \exists m_\varepsilon, \text{ s.t. } \forall m > m_\varepsilon \text{ and } n > m(\frac{1}{C_p} + \varepsilon),$
 $\exists f: \{0, 1\}^m \mapsto \{0, 1\}^n \text{ and } g: \{0, 1\}^n \mapsto \{0, 1\}^m, \text{ s.t. } \forall \mathbf{x} \in \{0, 1\}^m:$
$$\Pr_{z \leftarrow Z = (Z_1, \dots, Z_n)} [g(f(\mathbf{x}) \oplus z) \neq \mathbf{x}] \leq \varepsilon$$

for $Z_1, \dots, Z_n \text{ iid} \sim (1 - p, p)$.

- ▶ $C_p = 1 - h(p)$ — the **channel capacity**

$$p = .1 \implies C_p = 0.5310 > \frac{1}{2}$$

$$p = .25 \implies C_p \approx \frac{1}{5}$$

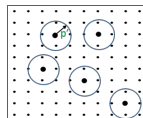
- ▶ Tight theorem
- ▶ We prove a weaker variant that holds w.h.p. over $\mathbf{x} \leftarrow \{0, 1\}^m$

Hamming distance

- ▶ For $\mathbf{y} = (y_1, \dots, y_n) \in \{0, 1\}^n$, let $|\mathbf{y}| = \sum_i y_i$ — Hamming weight of \mathbf{y}
- ▶ $|\mathbf{y} - \mathbf{y}'| = |\mathbf{y} \oplus \mathbf{y}'|$ — Hamming distance of \mathbf{y} from \mathbf{y}' ; # of places differ.

Proving the theorem

- ▶ Fix $p \in [0, \frac{1}{2}]$ and $\varepsilon > 0$, and let $m > m_\varepsilon$ and $n \geq m(\frac{1}{C_p} + \varepsilon)$, for m_ε to be determined by the analysis. (Recall $C_p = 1 - h(p)$).
- ▶ We show $\exists f: \{0, 1\}^m \mapsto \{0, 1\}^n$ and $g: \{0, 1\}^n \mapsto \{0, 1\}^m$, s.t.
 $\Pr_{\mathbf{x} \leftarrow \{0, 1\}^m} [g(f(\mathbf{x}) \oplus \mathbf{Z}) \neq \mathbf{x}] \leq \varepsilon$
- ▶ Main idea, find f s.t. $f(\mathbf{x})$ is $2p'n$ far from $f(\mathbf{x}')$ for all $\mathbf{x} \neq \mathbf{x}'$, for $p' > p$ to be determined later.



- ▶ Since $|\mathbf{Z}| \approx pn$, w.h.p. $|\mathbf{Z}| < p'$, and the noise won't hurt decoding.
- ▶ $g(y)$ returns $\operatorname{argmin}_{\mathbf{x}' \in \{0, 1\}^m} |y - f(\mathbf{x}')|$
- ▶ So it all boils down to finding f s.t.
 $\Pr_{\mathbf{x} \leftarrow \{0, 1\}^m; \mathbf{y} = f(\mathbf{x}) \oplus \mathbf{Z}} [|\mathbf{f}(\mathbf{x}) - \mathbf{y}| < \min_{\mathbf{x}' \in \{0, 1\}^m \setminus \{\mathbf{x}\}} |\mathbf{f}(\mathbf{x}') - \mathbf{y}|] \geq 1 - \varepsilon$
- ▶ We choose f uniformly at random (what does it mean?)
- ▶ Non-constructive proof
- ▶ Probabilistic method

Proving there exists good f

- ▶ For $y \in \{0, 1\}^n$, let $B_{p'}(y) = \{y' \in \{0, 1\}^n : |y' - y| \leq p'n\}$
- ▶ By weak law of large numbers, $\exists n' \in \mathbb{N}$ s.t. $\forall n \geq n'$ and $\forall \mathbf{x} \in \{0, 1\}^n$:
 $\Pr_{z \leftarrow Z} [(f(\mathbf{x}) \oplus z) \notin B_{p'}(f(\mathbf{x}))] \leq \frac{\varepsilon}{2}$ (for any fixed f)
- ▶ Fact (proved later): $b(p') = |B_{p'}(y)| \leq 2^{n \cdot h(p')}$

$$\Rightarrow \text{For all } \mathbf{x} \neq \mathbf{x}' \in \{0, 1\}^n: \Pr_f[f(\mathbf{x}') \in B_{p'}(f(\mathbf{x}))] = \frac{b(p')}{2^n} \leq \frac{2^{n \cdot h(p')}}{2^n} = 2^{-nC_{p'}}$$

$$\Rightarrow \forall \mathbf{x} \in \{0, 1\}^n: \Pr_f[\exists \mathbf{x}' \neq \mathbf{x} \in \{0, 1\}^m: f(\mathbf{x}') \in B_{p'}(f(\mathbf{x}))] \leq 2^{m-nC_{p'}}$$

$$\Rightarrow \exists f \text{ s.t. } \alpha_{m,n} = \Pr_{\mathbf{x} \leftarrow \{0,1\}^m} [\exists \mathbf{x}' \neq \mathbf{x} \in \{0, 1\}^m: f(\mathbf{x}') \in B_{p'}(f(\mathbf{x}))] \leq 2^{m-nC_{p'}}$$

$$\text{▶ } \alpha_{m,n} \leq \frac{\varepsilon}{2} \text{ for } n \geq \frac{m}{C_{p'}} - \log \varepsilon + 1 = m \left(\frac{1}{C_{p'}} + \frac{1 - \log \varepsilon}{m} \right)$$

$$\text{▶ Fix } p' > p \text{ such that } \frac{1}{C_{p'}} - \frac{1}{C_p} \leq \frac{\varepsilon}{2}$$

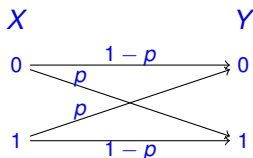
$$\Rightarrow \alpha_{m,n} \leq \frac{\varepsilon}{2} \text{ for } m \geq m' := \frac{2(1 - \log \varepsilon)}{\varepsilon} \text{ and } n \geq m \left(\frac{1}{C_p} + \frac{\varepsilon}{2} + \frac{1 - \log \varepsilon}{m} \right) \geq m \left(\frac{1}{C_p} + \varepsilon \right)$$

$$\text{▶ Hence, for } m > m_\varepsilon = \max\{m', n'\} \text{ and } n > m \left(\frac{1}{C_p} + \varepsilon \right), \text{ it holds that}$$

$$\Pr_{\mathbf{x} \leftarrow \{0,1\}^m} [g(f(\mathbf{x}) \oplus Z) \neq \mathbf{x}] \leq \varepsilon \quad \square$$

Why $C_p = 1 - h(p)$?

- ▶ Let $X \leftarrow \{0, 1\}$, $Z \sim (1 - p, p)$ and $Y = X \oplus Z$



- ▶ $I(X; Y) = H(Y) - H(Y|X) = H(Y) - H(Z) = 1 - h(p) = C_p$
- ▶ Received bit “gives” C_p information about transmitted bit
- ▶ Hence, to recover m bits, we need to send at least $m \cdot \frac{1}{C_p}$ bits

Size of bounding ball

Claim 2

For $p \in [0, \frac{1}{2}]$ and $n \in \mathbb{N}$: it holds that $\sum_{k=0}^{\lfloor pn \rfloor} \binom{n}{k} \leq 2^{n \cdot h(p)}$

Proof in a few slides (we already saw that $\binom{n}{pn} \approx 2^{n \cdot h(p)}$)

Corollary 3

For $y \in \{0, 1\}^n$ and $p \in [0, \frac{1}{2}]$, let $B_p(y) = \{y' \in \{0, 1\}^n : |y' - y| \leq pn\}$. Then $|B_p(y)| = \sum_{k=0}^{\lfloor pn \rfloor} \binom{n}{k} \leq 2^{n \cdot h(p)}$

Very useful estimation. Weaker variants follows by AEP or Stirling,

Tightness

- ▶ $X \leftarrow \{0, 1\}^m$, $Z = (Z_1, \dots, Z_n)$ where Z_1, \dots, Z_n iid $\sim (1 - p, p)$
- ▶ $\underbrace{X}_{m \text{ bits}} \rightarrow \underbrace{f(X)}_{n \text{ bits}} \rightarrow \underbrace{f(X) \oplus Z}_Y \rightarrow \underbrace{g(f(X) \oplus Z)}_{g(Y)}$
- ▶ Assuming $\Pr[g(Y) = X] \geq 1 - \varepsilon$, we show $nC_p \geq m(1 - \varepsilon) - 1$
- ▶ Compare to $nC_p > m(1 + \varepsilon C_p)$ in Thm 1
- ▶ Hence, $\lim_{\varepsilon \rightarrow 0} \frac{m}{n} = C_p$
- ▶ By Fano, $H(X|Y) \leq h(\varepsilon) + \varepsilon \log(2^m - 1) \leq 1 + \varepsilon m$
- ▶ $I(X; Y) = H(X) - H(X|Y) \geq m - \varepsilon m - 1 = m(1 - \varepsilon) - 1$
- ▶ $H(Y|X) = H(X, Y) - H(X) = H(X, Z) - H(X) = H(Z) = nh(p)$
- ▶ $I(X; Y) = H(Y) - H(Y|X) = n - nh(p)$
- ▶ Hence, $m(1 - \varepsilon) \leq I(X; Y) + 1 = n(1 - h(p)) + 1 = nC_p + 1$
- ▶ ...

General communication channel

$Q: [k] \mapsto [k]$ that channel (a probabilistic function)

$$p_{i,j} = \Pr[Q(i) = j]$$

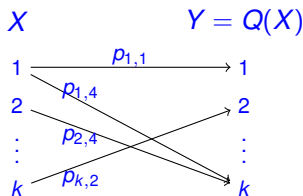
- ▶ $\mathbf{x} = (x_1, \dots, x_m) \in \{0, 1\}^m$
- ▶ Encoding function $f: \{0, 1\}^m \mapsto \{1, \dots, k\}^n$
- ▶ Decoding function $g: \{1, \dots, k\}^n \mapsto \{0, 1\}^m$
- ▶ $\mathbf{x} \xrightarrow{\text{encoding}} f(\mathbf{x}) \xrightarrow{\text{channel}} Q(f(\mathbf{x})) \xrightarrow{\text{decoding}} g(Q(f(\mathbf{x})))$
- ▶ We hope for $g(Q(f(\mathbf{x}))) = \mathbf{x}$

▶ **Channel capacity** $C_Q = \max_X I(X; Y)$

▶ The maximal information Y gives on X

▶ Shannon theorem: $\forall T$ and $\forall \varepsilon > 0$, $\exists m_\varepsilon$: $\forall m > m_\varepsilon$ and $\forall n > m(\frac{1}{C_Q} + \varepsilon)$:
 $\exists f, g$ as above s.t. $\Pr_Q[g(Q(f(\mathbf{x}))) \neq \mathbf{x}] \leq \varepsilon$, for all $\mathbf{x} \in \{0, 1\}^m$.

▶ Proof: similar lines to the binary case, but more subtle distribution for f



Discussion

- ▶ Tight result
- ▶ Non-constructive
- ▶ Coding theory: design **explicit** (and efficient) code achieving the same bounds
- ▶ Application: faulty communication, storage
- ▶ Combination of data compression and ECC

Part II

Combinatorial Applications

Movies

- ▶ 2^n people, $m = 3n$ movies.
- ▶ Every pair of movies was seen by at least 90% of the people
- ▶ Claim: there exist two people who saw exactly the same set of movies
- ▶ $X \leftarrow [2^n]$ — a random person
- ▶ $g_i(x) = \begin{cases} 1, & x \text{ saw movie } i; \\ 0, & \text{otherwise.} \end{cases}$
- ▶ $Y_i = g_i(X)$
- ▶ $\forall i, j: H(Y_i, Y_j) \leq H(.9, \frac{1}{3}, \frac{1}{3}, \frac{1}{3}) \leq \frac{2}{3}$
- ▶ $H(Y = (Y_1, \dots, Y_m)) \leq \sum_i^{m/2} H(Y_i, Y_{i+\frac{m}{2}}) < \frac{3n}{2} \cdot \frac{2}{3} = n = H(X)$
- ▶ Hence, X is not determined by Y

Why $H(X_1, \dots, X_n) \leq \sum_i H(X_i)$ so useful?

- ▶ $\mathcal{S} \subseteq \{0, 1\}^n$; $X = (X_1, \dots, X_n) \leftarrow \mathcal{S}$
- ▶ $\log |\mathcal{S}| = H(X) \leq \sum_i H(X_i)$ implies $|\mathcal{S}| \leq 2^{\sum_i H(X_i)}$
- ▶ If $\sum_i H(X_i)$ is small, then \mathcal{S} is small
 X_i are unbalanced, e.g., $\sim (.1, .9)$, implies $|\mathcal{S}| \leq 2^{n \cdot h(.1)} \leq 2^{n/2}$
- ▶ \mathcal{S} is large implies $\sum_i H(X_i)$ is large, hence most X_i are almost balanced
- ▶ $|\mathcal{S}| \geq \frac{1}{2} 2^n$ implies $\mathbb{E}_{i \leftarrow [n]} [H(X_i)] \geq 1 - \frac{1}{n}$
- ▶ Most X_i are close to uniform

Hamming ball

- ▶ $p \leq \frac{1}{2}$; $\mathcal{S} = \{(a_1, \dots, a_n) \in \{0, 1\}^n : \sum_i a_i \leq pn\}$
- ▶ $|\mathcal{S}| = \sum_{k=0}^{\lfloor pn \rfloor} \binom{n}{k}$
- ▶ $X = (X_1, \dots, X_n) \leftarrow \mathcal{S}$
- ▶ $\sum_i X_i \leq pn \implies \mathbb{E}[\sum X_i] \leq pn$, and by symmetry $\mathbb{E}[X_i] \leq p$ for every i
- ▶ $\forall i, j: \Pr[X_i = 1] = \Pr[X_j = 1] \leq p$

$\implies H(X_i) \leq h(p)$ for every i

$\implies |\mathcal{S}| \leq 2^{nh(p)}$

$\implies \sum_{k=0}^{\lfloor pn \rfloor} \binom{n}{k} \leq 2^{nh(p)}$

▶ ...

Application

- ▶ X_1, \dots, X_n iid uniform bits (i.e., $\sim (\frac{1}{2}, \frac{1}{2})$)
- ▶ $\Pr[\sum_i X_i \leq pn] = \Pr[(X_1, \dots, X_n) \in \mathcal{S}] \leq 2^{nh(p)} \cdot 2^{-n} = 2^{-n(1-h(p))}$
- ▶ Very useful inequality. No Chernoff, just IT

Isoperimetric inequality

- ▶ $\mathcal{S} \subseteq \{0, 1\}^n$
- ▶ Edges of \mathcal{S} — $E = \{(u, v) \in \mathcal{S} : |u - v| = 1\}$

Theorem 4

$$|E| \leq \frac{1}{2} \cdot |\mathcal{S}| \cdot \log |\mathcal{S}|$$

- ▶ Equality if \mathcal{S} is “face” : $\mathcal{S} = \{(\mathbf{x}, \mathbf{y}) : \mathbf{y} \in \{0, 1\}^d\}$ for some $\mathbf{x} \in \{0, 1\}^{n-d}$
- ▶ Example: \mathcal{S} is a **face** of the 3-dimensional cube
 $n = 3, |\mathcal{S}| = 4$, implies $|E| \leq \frac{1}{2} \cdot 4 \cdot \log 4 = 4$
- ▶ E_i — edges of E in **direction** i ($E = \bigsqcup_{i \in [n]} E_i$)
- ▶ $X = (X_1, \dots, X_n) \leftarrow \mathcal{S}$ and $X_{-i} = (X_1, X_2, \dots, X_{i-1}, X_{i+1}, \dots, X_n)$

Lemma 5

$$H(X_i | X_{-i}) = \frac{2|E_i|}{|\mathcal{S}|}$$

Proving **Thm 4**:

$$\begin{aligned} \log |\mathcal{S}| &= H(X_1, \dots, X_n) = H(X_1) + H(X_2 | X_1) + \dots + H(X_n | X_1, X_2, \dots, X_{n-1}) \\ &\geq H(X_1 | X_{-1}) + H(X_2 | X_{-2}) + \dots + H(X_n | X_{-n}) = \sum_i \frac{2|E_i|}{|\mathcal{S}|} = \frac{2|E|}{|\mathcal{S}|}. \quad \square \end{aligned}$$

Proving Lemma 5

We prove for $i = 1$

- ▶ $\mathcal{S} \subseteq \{0, 1\}^n$; $X = (X_1, \dots, X_n) \leftarrow \mathcal{S}$
- ▶ $E = \{(u, v) \in \mathcal{S} : |u - v| = 1\}$ and E_1 contains edges of E in direction 1
- ▶ $\mathcal{S}_{-1} = \{\mathbf{y} \in \{0, 1\}^{n-1} : \exists x \in \{0, 1\} \text{ s.t. } (x, \mathbf{y}) \in \mathcal{S}\}$.
(\mathcal{S} projected on $(2, \dots, n)$)
- ▶ $\mathcal{S}_{-1}^e = \{\mathbf{y} \in \{0, 1\}^{n-1} : (0, \mathbf{y}), (1, \mathbf{y}) \in \mathcal{S}\}$ and $\mathcal{S}_{-1}^{\neg e} = \mathcal{S}_{-1} \setminus \mathcal{S}_{-1}^e$
- ▶ $|\mathcal{S}| = 2|\mathcal{S}_{-1}^e| + |\mathcal{S}_{-1}^{\neg e}|$
- ▶ $|E_1| = |\mathcal{S}_{-1}^e|$
- ▶ $H(X|X_{-1}) = \Pr[X_{-1} \in \mathcal{S}_{-1}^e] \cdot 1 = \frac{2|\mathcal{S}_{-1}^e|}{2|\mathcal{S}_{-1}^e| + |\mathcal{S}_{-1}^{\neg e}|} = \frac{2|E_1|}{|\mathcal{S}|}$ □
- ▶ ...